Query Rewriting with Symmetric Constraints

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Abstract. We address the problem of answering queries using expressive *symmetric* inter-schema constraints which allow to establish mappings between several heterogeneous information systems. This problem is of high relevance to data integration, as symmetric constraints are essential for dealing with true concept mismatch and are generalizations of the kinds of mappings supported by both local-as-view and global-as-view approaches that were previously studied in the literature. Moreover, the flexibility gained by using such constraints for data integration is essential for virtual enterprise and e-commerce applications. We first discuss resolution-based methods for computing maximally contained rewritings and characterize computability aspects. Then we propose an alternative but semantically equivalent perspective based on a generalization of results relating to the database-theoretic problem of answering queries using views. This leads to a fast query rewriting algorithm, which has been implemented and experimentally evaluated.

1 Introduction

This paper addresses the *query rewriting problem* in semantic data integration in a very general form, as a proper generalization of the well-known local-as-view (e.g., [17,11,2]) and global-as-view approaches (e.g., [10,1,5]). To start somewhere, we focus on the relational case. Given a conjunctive query Q, we attempt to find a *maximally contained rewriting* in terms of a set of distinguished *source predicates* $\mathbf{S}$ only – under a given set of constraints, the positive relational queries as the output query language (i.e., a rewriting is a set of conjunctive queries), and the classical logical semantics. Under the classical semantics,

1. for each conjunctive query Q' in the maximally contained rewriting of Q, the constraints taken as a logical theory imply $Q \supseteq Q'$ and Q' uses predicates of $\mathbf{S}$ only, and
2. for each conjunctive query Q'' over predicates in $\mathbf{S}$ only for which the constraints imply that $Q \supseteq Q''$, there is a conjunctive query Q' in the rewriting such that the constraints imply $Q' \supseteq Q''$.

We support inter-schema constraints in the form of what we call *Conjunctive Inclusion Dependencies* (cind's), containment relationships between conjunctive queries. We refer to cind's as *symmetric* because they are syntactically symmetric with respect to the inclusion resp. implication symbol, while for instance materialized view definitions used for local-as-view query rewriting are not.

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Example 1.1. Consider a conjunctive query

$$Q(x_1) \leftarrow \text{parent}(x_1, x_2),\ \text{parent}(x_2, x_3),\ \text{parent}(x_3, x_4).$$

asking for great-grandparents in terms of a schema which contains a predicate "parent". Now let "parent" be a "logical" relation to which database relations first need to be mapped before any queries over it can be answered. Let "parent" be conceived to represent parent-child relationships between living persons only. We want to rewrite Q into a query over source predicates in $\mathbf{S} = \{\text{grandparent}, \text{alive}\}$, where "grandparent" contains grandparent relationships between persons of which the grandparents may possibly have deceased and the source relation "alive" holds persons still alive. We may assert the cind

$$\{\langle x, z\rangle \mid \exists y : \text{parent}(x, y) \wedge \text{parent}(y, z)\} \supseteq \{\langle x, z\rangle \mid \text{grandparent}(x, z) \wedge \text{alive}(x)\}$$

between the schema of "parent" and $\mathbf{S}$, i.e., a mapping formulated as a containment relationship between conjunctive queries. Then,

$$Q'(x_1) \leftarrow \text{grandparent}(x_1, x_3),\ \text{alive}(x_1),\ \text{grandparent}(x_3, z),\ \text{alive}(x_3).$$

is the maximally contained rewriting of Q, i.e. the largest query logically contained in Q with respect to the constraint given (and thus cannot return any wrong answers) that can be computed by only considering query and constraint, but not the data in the relations of $\mathbf{S}$, and which only uses relations from $\mathbf{S}$. □

Note that the problem we attack in this paper is different from work on query answering using integrity constraints [8,9,12] or the problem of optimizing queries in mediated systems. Our constraints are meant to encode mappings between the schemata of several information systems as in [4,3]; as such, they compare to materialized view *definitions* in the local-as-view approach rather than to classical *integrity constraints*.

cind's may represent mappings between complex networks of information systems in which each information system may have its own schema and make use of integrated data. We are thus not restricted to a single layer of constraints (which map sources against a so-called "global" schema) as in the local-as-view approach. Starting from a set of actual source relations which contain data, views or "logical relations" can be defined and used to extend any of the schemata. To facilitate practical usability, a logical relation may map to sources through several indirections. Such an architecture is essential for large and open environments such as experienced in the modern e-economy [4,14].

The relevance of query rewriting with symmetric constraints stems from the observation that both local-as-view and global-as-view approaches to data integration are unable to deal with concept mismatch requiring mappings between pairs of complex query expressions. This problem is particularly common in complex technical domains [14].

Example 1.2. Assume that we have a number of sources holding information about computer models as-built (in fact, we only consider computer mainboards

in this short example), their constituent parts, and technical specifications. We want to integrate these sources against a reference design view **R** with predicates "mb" (for mainboards), "cpu", "cache", "conn" (for connections or part-of relationships between component parts), and possibly others. Let a source schema **S** represent mainboards using a CPU model called "p1" which has an on-chip cache. We cannot directly map cache components in this example, but we can map characteristics represented by both schemata (say, MHz rates r and cache sizes s) to mainboards and CPUs and mediate useful queries over **R**. We encode the desired mapping using the cind

$$\{\langle x,y,r,s\rangle \mid \exists z : \mathbf{R}.\mathrm{mb}(x),\ \mathbf{R}.\mathrm{conn}(x,y),\ \mathbf{R}.\mathrm{conn}(x,z),\ \mathbf{R}.\mathrm{cpu}(y,r),\ \mathbf{R}.\mathrm{cache}(z,s)\} \supseteq \{\langle x,y,r,s\rangle \mid \mathbf{S}.\mathrm{mb}(x),\ \mathbf{S}.\mathrm{conn}(x,y),\ \mathbf{S}.\mathrm{p1}(y,r,s)\}$$ □

Local-as-view and global-as-view approaches assume that mediated schemata can be designed beforehand using intuitions concerning the likely sources to be added to an integration system later. Both large data integration settings and changing requirements render such an assumption unsustainable (see [14]). Our approach allows to "patch" local-as-view or global-as-view integration systems when sources need to be integrated whose particularities have not been foreseen when designing the schemata against which data are to be integrated.

The only previous work dealing with symmetric constraints is the description logics approach to data integration (e.g., [3,4]), which, however, requires high-complexity reasoning over the data (thus, there is a scalability issue) and to import all data to be integrated into the description logics reasoner. This is usually not an option in open e-economy or WWW data integration environments. Solving the integration problem on the level of queries and mappings *only* is essential for being able to deal with large amounts of data and restricted (e.g., screen-scraping) query interfaces.

Contributions and Structure. After some preliminaries (Section 2), we discuss a simple resolution-based method for generating rewritings and provide characterizations of the main theoretical properties of our problem in Section 3. Unfortunately, positive (and thus non-recursive) maximally contained rewritings may be infinite and the major decision problems (such as the non-emptiness or boundedness of the result) are undecidable. However, given that the predicate dependency graph (with respect to the inclusion direction) of a set of constraints is acyclic, we can guarantee to find the maximally contained rewritings, which are finite. The acyclic case is a proper generalization of both local-as-view and global-as-view approaches.

In Section 4, we propose an alternative algorithm for computing maximally contained rewritings which is based on a generalization of the MiniCon Algorithm [18] for the problem of answering queries using views, and demonstrate its soundness and completeness. When using this algorithm, all intermediate rewritings are guaranteed to be function-free and thus conjunctive queries. Because of that, one can make use of classical database techniques for optimizing the rewriting process. Section 5 presents refinements of the algorithm of Section 4, which we have implemented in a practical system.

We evaluate our implementation, which is publicly available, experimentally in Section 6. It turns out that it scales to thousands of constraints and realistic applications. Section 7 concludes with a discussion of our new algorithm.

2 Preliminaries

We define a *conjunctive inclusion dependency* (cind) as a constraint of the form $Q_1 \subseteq Q_2$ where Q_1 and Q_2 are conjunctive queries of the form

$$\{\langle x_1, \ldots, x_n\rangle \mid \exists x_{n+1} \ldots x_m : (p_1(\bar{X}_1) \wedge \ldots \wedge p_k(\bar{X}_k))\}$$

with a set of *distinct*[1] unbound variables $x_1, \ldots, x_n$, without arithmetic comparisons, but possibly with constants. We may write $\{Q_1 \equiv Q_2\}$ as a short form of $\{Q_1 \subseteq Q_2,\ Q_1 \supseteq Q_2\}$. The *normal form* $NF(\Sigma)$ of a set Σ of cind's – i.e., Σ taken as a logical formula transformed into (implication) normal form – is a set of Horn clauses of a simple pattern. Every cind σ of the form $Q_1 \subseteq Q_2$ with

$$Q_1 = \{\langle x_1, \ldots, x_n\rangle \mid \exists x_{n+1} \ldots x_m : v_1(\bar{X}_1) \wedge \ldots \wedge v_k(\bar{X}_k)\}$$
$$Q_2 = \{\langle y_1, \ldots, y_n\rangle \mid \exists y_{n+1} \ldots y_{m'} : p_1(\bar{Y}_1) \wedge \ldots \wedge p_{k'}(\bar{Y}_{k'})\}$$

translates to k' Horn clauses $p_i(\bar{Z}_i) \leftarrow v_1(\bar{X}_1) \wedge \ldots \wedge v_k(\bar{X}_k))$. where each $z_{i,j}$ of $\bar{Z}_i$ is determined as follows: If $z_{i,j}$ is a variable y_h with $1 \leq h \leq n$, replace it with x_h. If $z_{i,j}$ is a variable y_h with $n < h \leq m'$, replace it with Skolem function $f_{\sigma,y_h}(x_1, \ldots, x_n)$ (the subscript assures that the Skolem functions are unique for a given constraint and variable).

Example 2.1. Let σ be the cind

$$\{\langle y_1, y_2\rangle \mid \exists y_3 : p_1(y_1, y_3) \wedge p_2(y_3, y_2)\} \supseteq \{\langle x_1, x_2\rangle \mid \exists x_3 : v_1(x_1, x_2) \wedge v_2(x_1, x_3)\}$$

Then, $NF(\{\sigma\})$ is

$$p_1(x_1, f_{\sigma,y_3}(x_1, x_2)) \leftarrow v_1(x_1, x_2) \wedge v_2(x_1, x_3).$$
$$p_2(f_{\sigma,y_3}(x_1, x_2), x_2) \leftarrow v_1(x_1, x_2) \wedge v_2(x_1, x_3).$$ □

Whenever a cind translates into a function-free clause in normal form, we write it in datalog notation. This is the case for cind's $\{\langle \bar{X}\rangle \mid p(\bar{X})\} \supseteq Q$, i.e., where the subsumer queries are $\exists$-free single-literal queries.

The dependency graph of a set C of Horn clauses is the directed graph constructed by taking the predicates of C as nodes and adding, for each clause in C, an edge from each of the body predicates to the head predicate. The diameter of a directed acyclic graph is the longest directed path occurring in it. The dependency graph of a set of cind's Σ is the dependency graph of the logic program $NF(\Sigma)$. A set of cind's is *cyclic* if its dependency graph is cyclic. An acyclic set Σ of cind's is called *layered* if the predicates appearing in Σ can be partitioned into n disjoint sets $P_1, \ldots, P_n$ such that there is an index i for each cind $\sigma : Q_1 \subseteq Q_2 \in \Sigma$ such that $Preds(Body(Q_1)) \subseteq P_i$ and $Preds(Body(Q_2)) \subseteq P_{i+1}$ and $Sources = P_1$.

[1] Note that if we would not require unbound variables in constituent queries to be distinct, the transformation into normal form would result in Horn clauses with equality atoms as heads.

3 Query Containment and Rewriting

Let us begin with a straightforward remark on the containment problem for conjunctive queries under a set of cind's Σ, which, since they are themselves containment relationships between conjunctive queries, is the implication problem for this type of constraint. If we want to check a containment $\{\langle \bar{X} \rangle \mid \exists \bar{Y} : \phi(\bar{X}, \bar{Y})\} \supseteq \{\langle \bar{X} \rangle \mid \exists \bar{Z} : \psi(\bar{X}, \bar{Z})\}$ of two conjunctive queries under Σ by refutation (without loss of generality, we assume $\bar{Y}$ and $\bar{Z}$ to be disjoint and the unbound variables in the two queries above to be the same, $\bar{X}$), we have to show $\Sigma, \neg(\forall \bar{X} : (\exists \bar{Y} : \phi(\bar{X}, \bar{Y})) \leftarrow (\exists \bar{Z} : \psi(\bar{X}, \bar{Z}))) \vDash \bot$ i.e. the inconsistency of the constraints and the negation of the containment taken together. In normal form, ψ becomes a set of ground facts where all variables have been replaced *one-to-one* by new constants and ϕ becomes a clause with an empty head, where all distinguished variables x_i have been replaced by constants also used in ψ.

Example 3.1. For proving

$$\Sigma \vDash \{\langle x_1, x_2 \rangle \mid \exists x_3 : (p_1(x_1, x_3) \wedge p_2(x_3, x_2))\} \supseteq \{\langle y_1, y_2 \rangle \mid \exists y_3 : (r_1(y_1, y_3) \wedge r_2(y_3, y_2))\}$$

for a set of cind's Σ, we have to create the logic program

$$\mathcal{P} := NF(\Sigma) \cup \{ \leftarrow p_1(\alpha_1, x_3) \wedge p_2(x_3, \alpha_2). \quad r_1(\alpha_1, \alpha_3) \leftarrow . \quad r_2(\alpha_3, \alpha_2) \leftarrow . \}$$

where $\alpha_1, \alpha_2, \alpha_3$ are constants not appearing elsewhere. By the correctness of resolution for logic programs, the containment above holds iff there is a refutation of the goal $\leftarrow p_1(\alpha_1, x_3) \wedge p_2(x_3, \alpha_2).$ with the remaining clauses in $\mathcal{P}$. □

Definition 3.1. A set of conjunctive queries $\mathcal{Q}$ is a maximally contained positive rewriting of a conjunctive query Q with respect to a set of cind's Σ and a set of source predicates $\mathbf{S}$ iff

1. for each $Q' \in \mathcal{Q}$, $\Sigma \vDash Q \supseteq Q'$ and $Preds(Q') \subseteq \mathbf{S}$ and
2. for each conjunctive query Q'' with $\Sigma \vDash Q \supseteq Q''$ and $Preds(Q'') \subseteq \mathbf{S}$, there is a $Q' \in \mathcal{Q}$ such that $\Sigma \vDash Q \supseteq Q' \supseteq Q''$. □

In the finite case, a minimal[2] such set $\mathcal{Q}$ is of course unique up to reordering and variable renaming.

For simplicity, we assume that no source predicates appear in any heads of Horn clauses in $NF(\Sigma)$ throughout this paper. This does not cause any loss of generality, since we can always replace a source predicate that violates this assumption by a new virtual predicate in all cind's and then add a cind that maps the source predicate to that new virtual predicate.

Informally, we can obtain such a maximally contained rewriting by a method based on SLD resolution in the following way. Given a conjunctive query Q, a set of cind's Σ, and a set of source predicates $\mathbf{S}$, we first create a logic program

[2] A set of conjunctive queries is minimal if and only if the constituent queries are individually minimal and pairwise non-redundant.

$NF(\Sigma)$ and add a unit clause $s(\bar{X}) \leftarrow .$ (with a tuple $\bar{X}$ of *distinct* variables) for each predicate $s \in \mathbf{S}$. Then we try to refute the body of Q. (Differently from what we do for containment, we do not freeze any variables.) If we have found a refutation with a most general unifier θ, we collect the unit clauses used and create a Horn clause with $\theta(Head(Q))$ as head and the application of θ to the copies of unit clauses involved in the proof as body. If this clause is function-free, we output it. After that *we go on* as if we had not found a "proof" to compute more rewritings. Given an appropriate selection rule or a breath-first strategy for computing derivations, it is easy to see that this method will compute a maximally contained rewriting of Q in terms of multi-sets of conjunctive queries in the sense that for each conjunctive query Q'' contained in Q, a subsumer Q' will eventually be produced s.t. $\Sigma \models Q \supseteq Q' \supseteq Q''$. See Example 4.2 for query rewriting by an altered refutation proof.

Computability and Complexity.

Theorem 3.1. Let Σ be a set of cind's, $\mathbf{S}$ a set of predicates, and Q and Q' be conjunctive queries. Then the following problems are undecidable:

1. $\Sigma \models Q \subseteq Q'$, the containment problem.
2. $\exists Q' : \Sigma \models Q \supseteq Q'$ s.t. $Preds(Q') \subseteq \mathbf{S}$, i.e. it is undecidable whether the maximally contained rewriting of a conjunctive query Q w.r.t. Σ and $\mathbf{S}$ is nonempty (that is, it contains at least one conjunctive query). □

Moreover, the boundedness problem for maximally contained positive rewritings is undecidable, as any datalog program can be written as a set of cind's.

We next give an intuition for the undecidability results of Theorem 3.1. Post's Correspondence Problem (PCP, see e.g. [20]), a simple and well-known undecidable problem, is defined as follows. Given nonempty words $x_1, \ldots, x_n$ and $y_1, \ldots, y_n$ over the alphabet $\{0, 1\}$, the problem is to decide whether there are indexes $i_1, \ldots, i_k$ (with $k > 0$) s.t. $x_{i_1} x_{i_2} \ldots x_{i_k} = y_{i_1} y_{i_2} \ldots y_{i_k}$. Pairs of words $\langle x_i, y_i \rangle$ are also called *dominos*. In the following example, we show, by an example, an encoding of PCP in terms of our query rewriting problem.

Example 3.2. Let s be a source, $q \leftarrow inc(0,0).$ a boolean query, and

$$\mathrm{inc}(x,y) \leftarrow \mathrm{one}(x,x_1),\ \mathrm{zero}(x_1,x_2),\ \mathrm{one}(x_2,x_3),\ \mathrm{one}(y,y_1),\ \mathrm{inc}(x_3,y_1). \quad (1)$$

$$\mathrm{inc}(x,y) \leftarrow \mathrm{one}(x,x_1),\ \mathrm{zero}(y,y_1),\ \mathrm{one}(y_1,y_2),\ \mathrm{one}(y_2,y_3),\ \mathrm{one}(y_3,y_4),\ \mathrm{zero}(y_4,y_5),\ \mathrm{inc}(x_1,y_5). \quad (2)$$

$$\mathrm{inc}(x,y) \leftarrow \mathrm{dec}(x,y). \quad (3)$$

$$\{\langle x,y\rangle \mid \mathrm{dec}(x,y)\} \subseteq \{\langle x,y\rangle \mid \exists x_1,y_1 : \mathrm{zero}(x,x_1) \wedge \mathrm{zero}(y,y_1) \wedge \mathrm{dec}(x_1,y_1)\} \quad (4)$$

$$\{\langle x,y\rangle \mid \mathrm{dec}(x,y)\} \subseteq \{\langle x,y\rangle \mid \exists x_1,y_1 : \mathrm{one}(x,x_1) \wedge \mathrm{one}(y,y_1) \wedge \mathrm{dec}(x_1,y_1)\} \quad (5)$$

$$\mathrm{dec}(0,0) \leftarrow s. \quad (6)$$

six cind's of which the leading two stand for the instance

$$\mathcal{I} = \{\langle x_1 = 101,\ y_1 = 1\rangle, \langle x_2 = 1,\ y_2 = 01110\rangle\}$$

and the remaining four constitute the core PCP encoding. The constraints (1) and (2) "guess" two words represented as chains of "one" and "zero" atoms by the nondeterminism by which resolution (or MCD rewriting using Algorithm 4.1, for that matter) chooses a clause to resolve an "inc" atom, (3) finalizes the guess phase, constraints (4) and (5) "check" whether the two words are equal (which indicates the existence of a solution to the PCP problem) by proceeding from the right to the left, and constraint (6) "terminates" if the search was successful.

For showing the PCP instance $\mathcal{I}$ satisfiable, one can compute a contained rewriting by applying the constraints in the following order (we only describe the proof but no dead-end branches): (guess phase) (1), (2), (1), (3), (check phase) (5), (4), (5), (5), (5), (4), (5), (termination) (6)[3]. We find a solution $x_1x_2x_1 = y_1y_2y_1 = 1011101$ to $\mathcal{I}$. Generally, a PCP instance is satisfiable iff the maximally contained rewriting is $\{q \leftarrow s.\}$. (Furthermore, a PCP instance is satisfiable iff $\Sigma \vDash \{\langle\rangle \mid inc(0,0)\} \supseteq \{\langle\rangle \mid s\}$.) □

For the important case that Σ is acyclic, the above problems are decidable (and those of Theorem 3.1 are *NEXPTIME*-complete).

Theorem 3.2. Let Σ be an acyclic set of cind's and Q and Q' be conjunctive queries. Then the containment problem $\Sigma \vDash Q \subseteq Q'$ and the query rewriting problem for conjunctive queries (under acyclic sets of cind's) are NEXPTIME-complete. □

Membership in NEXPTIME follows from the more general result on nonrecursive logic programming [7,21] and hardness can be shown by a modification of the reduction from the TILING problem of [7]. For lack of space, we have refer to [14] for a proof.

4 Generalizing Local-as-View Rewriting

The results of this section generalize from work on algorithms for the problem of answering queries using views [16], for instance the Bucket Algorithm [17], the Inverse Rules Algorithm [8], OCCAM [15], the Unification-join Algorithm [19], and particularly the MiniCon Algorithm [18]. For space reasons, we introduce necessary notions as needed and refer to [18] for a discussion and comparison of such algorithms.

We adapt the notion of MiniCon descriptions [18] to our framework based on query rewriting with cind's decomposed into Horn clauses.

Definition 4.1. (Inverse MiniCon Description). Let Q be a conjunctive query with $n = |Body(Q)|$ and Σ be a set of cind's. An (inverse) MiniCon description for Q is a pair of a tuple $\langle c_1, \ldots, c_n\rangle \in (NF(\Sigma) \cup \{\epsilon\})^n$ with at least one $c_i \neq \epsilon$ and a substitution θ that satisfies the following two conditions.

[3] One can easily verify this proof using the intuition of fully replacing parts of (intermediate) goals by subsumed queries of cind's whose subsumer queries fully match those parts. Due to the special structure of the cind's, at any point, all MCDs are "isomorphic" to some subsumer query of a cind.

1. For the most general unifier $\theta \neq fail$ arrived at by unifying the heads of all the $c_i \neq \epsilon$ with $Body_i(Q)$, the unfolding of Q and $\langle c_1, \ldots, c_n \rangle$ under θ is function-free and
2. there is no tuple $\langle c'_1, \ldots, c'_n \rangle \in \{c_1, \epsilon\} \times \ldots \times \{c_n, \epsilon\}$ with fewer entries different from ϵ than in $\langle c_1, \ldots, c_n \rangle$, such that the unfolding of Q with $\langle c'_1, \ldots, c'_n \rangle$ is function free. □

Example 4.1. Consider again the query Q and the constraint (which we now call σ) of Example 1.1. $NF(\{\sigma\})$ is

$$c_1 : \ \text{parent}(x, f_{\sigma,y}(x,z)) \leftarrow \text{grandparent}(x,z) \wedge \text{alive}(x).$$
$$c_2 : \ \text{parent}(f_{\sigma,y}(x,z), z) \leftarrow \text{grandparent}(x,z) \wedge \text{alive}(x).$$

We have two MCDs, $\langle\langle c_1, c_2, \epsilon\rangle, \theta\rangle$ with the unifier

$$\theta = \{[x_1/x^{(1)}],\ [x_2/f_{\sigma,y}(x^{(1)}, z^{(1)})],\ [x_2/f_{\sigma,y}(x^{(2)}, z^{(2)})],\ [x_3/z^{(2)}]\}$$

and $\langle\langle \epsilon, \epsilon, c_1\rangle, \theta'\rangle$ with $\theta' = \{[x_3/x^{(3)}], [x_4/f_{\sigma,y}(x^{(3)}, z^{(3)})]\}$. Note that $\langle c_1, c_2, c_1\rangle$ violates condition 2 of Definition 4.1, while all other MCD candidates violate condition 1. □

Note that the inverse MiniCon descriptions of Definition 4.1 exactly coincide with the MCDs of [18] in the local-as-view case. Algorithm 4.1 shown below can easily be reformulated so as to use a slight generalization of the notation of [18] to cover clause bodies consisting of several atoms. That way, one can even escape the need to transform cind's into Horn clauses and can reason completely without the introduction of function terms. However, to support the presentation of our results (particularly the equivalence proof of the following section), we do not follow this path in this paper.

Algorithm 4.1 (Query rewriting with MCDs).
Input. A conjunctive query Q, a set of cind's Σ, and a set **S** of source predicates
Output. A maximally contained rewriting of Q w.r.t. Σ and **S**

$Qs := [Q]$;
while Qs is not empty **do** {
  $[Q, Qs] := Qs$;
  if $Preds(Q) \subseteq$ **S then output** Q;
  else {
    $M :=$ compute the set of all inverse MCDs for Q and Σ;
    for each $\langle\langle c_1, \ldots, c_n\rangle, \theta\rangle \in M$ **do** {
      $Q' :=$ unfold$(Q, \theta, \langle c_1, \ldots, c_n\rangle)$;
      $Qs := [Qs, Q']$;
} } } □

In Algorithm 4.1, maximally contained rewritings of a conjunctive query Q are computed by iteratively unfolding queries with *single* MiniCon descriptions[4]

[4] In this respect, Algorithm 4.1 differs from the MiniCon algorithm for the problem of answering queries using views, where MCDs are packed so as to rewrite all body atom at once.

until a rewriting contains only source predicates in its body. In order to handle cyclic sets of cind's (and attain completeness), we manage intermediate rewritings using a queue and, consequently, follow a breath-first strategy.

The function "unfold" accepts a conjunctive query Q with $|Body(Q)| = n$, a unifier θ and a tuple of n Horn clauses or ϵ s.t. if $c_i \neq \epsilon$, θ unifies $Body_i(Q)$ with $Head(c_i)$. It produces a new clause from Q (which in fact is again guaranteed to be function-free and thus a conjunctive query) by replacing $Head(Q)$ by $\theta(Head(Q))$ and each of the non-source body atoms $Body_i(Q)$, with $c_i \neq \epsilon$, by $\theta(Body(c_i))$. (i.e. after applying substitutions from the unifier). If $c_i = \epsilon$, $Body_i(Q)$ is replaced by $\theta(Body_i(Q))$. Of course, for each MCD $\langle\langle \ldots, c_i, \ldots, c_j, \ldots\rangle, \theta\rangle$ we have $\theta(Body(c_i)) = \theta(Body(c_j))$, and thus only one rule body needs to be added for each MCD during unfolding.

Theorem 4.2. Let Q be a conjunctive query, Σ be a set of cind's, and $\mathbf{S}$ be a set of "source" predicates. Then, for each conjunctive query Q' with $Preds(Q') \subseteq \mathbf{S}$ we have $\Sigma \models Q \supseteq Q'$ iff Algorithm 4.1 eventually computes a conjunctive query Q'' with $Preds(Q'') \subseteq \mathbf{S}$ and $\Sigma \models Q \supseteq Q'' \supseteq Q'$. □

In other words, Algorithm 4.1 enumerates the maximally contained positive rewriting of Q under Σ in terms of $\mathbf{S}$. Note that given our requirement that predicates of $\mathbf{S}$ do not appear in heads of $NF(\Sigma)$, we of course have $\Sigma \models Q'' \supseteq Q'$ iff $Q'' \supseteq Q'$ (that is, classical conjunctive query containment [6]).

It is easy to see that the rewriting process of Algorithm 4.1 simply is equivalent to resolution where only some of the subgoals of a goal may be rewritten in a single step and each intermediate rewriting has to be function-free. Every proof generated by Algorithm 4.1 is thus a correct resolution proof. Thus,

Lemma 4.1. (Soundness of Algorithm 4.1) Let Q be a conjunctive query, Σ a set of cind's, and $\mathbf{S}$ a set of source predicates. Then, for each conjunctive query Q' generated by Algorithm 4.1 for Q, Σ, and $\mathbf{S}$, we have $\Sigma \models Q \supseteq Q'$ and $Preds(Q') \subseteq \mathbf{S}$. □

Completeness is a consequence of the following result.

Lemma 4.2. Let $\mathcal{P}$ be a resolution proof establishing a logically contained rewriting of a conjunctive query Q under a set of cind's Σ. Then, there is always a proof $\mathcal{P}'$ establishing the same contained rewriting such that each intermediate rewriting is function-free. □

Proof. Let us assume that each new subgoal a derived using resolution receives an identifying index $idx(a)$. Then, given the proof $\mathcal{P}$, there is a unique *next premise to be applied* $c_{idx(a)}$ out of the Horn clauses in $NF(\Sigma)$ for each subgoal a. This is the Horn clause from our constraints base that will be unfolded with a to resolve it in $\mathcal{P}$.

Note that the proof $\mathcal{P}$ is fully described by some unique indexing $idx(a)$ of all subgoals a appearing in the proof (while we do not need to know or remember the atoms themselves), the clauses $c_{idx(a)}$, and a specification of which indexes the subgoals in the bodies of these clauses are attributed with when they are unfolded with subgoals.

In our original proof $\mathcal{P}$, each subgoal a of a goal is rewritten with $c_{idx(a)}$ in each step, transforming g_0, the body of Q and at the same time the initial goal, via $g_1, \ldots, g_{n-1}$ to g_n, the body of the resulting rewriting. We maintain the head of Q separately across resolution steps and require that variables in the head are not unified with function terms, but apply other unifications effected on the variables in the goals in parallel with the rewriting process. Already $\mathcal{P}$ must assure *at any step* that no variable from the head of Q is unified with a function term, as otherwise no conjunctive query can result.

We know that resolution remains correct no matter in which order the next due resolution steps $c_{idx(a)}$ are applied to the subgoals, and that we even may unfold, given e.g. a goal with two atoms, the first goal and then a subgoal from the unfolding of that first goal (and may do that any finite number of times) before we unfold our second original subgoal.

Coming back to deriving a function-free proof starting from $\mathcal{P}$, all we now have to show is that at any intermediate step of a resolution proof with cind's, a nonempty set of subgoals $X = \{a_{i_1}, \ldots, a_{i_k}\} \subseteq g_i$ of the function-free intermediate goal g_i exists such that, when only these subgoals are unfolded with their next due premises to be applied $c_{idx(a_{i_1})}, \ldots, c_{idx(a_{i_k})}$, the overall new goal g_{i+1} produced will be function-free[5]. The emphasis here lies on finding a *nonempty* such set X, as the empty set automatically satisfies this condition. If we can guarantee that such a nonempty set always exists until the function-free proof has been completed, our lemma is shown.

Let there be a dependency graph $G_{g_i} = \langle V, E \rangle$ for each (intermediate) goal g_i with the subgoals as vertices and a directed edge $\langle a, b \rangle \in E$ iff a contains a variable v that is unified with a function term $f(\bar{X})$ in $Head(c_{idx(a)})$ and v appears in b and is unified with a variable (rather than a function term with the same function symbol) in $Head(c_{idx(b)})$. (Intuitively, if there is an edge $\langle a, b \rangle \in E$, then b must be resolved *before* a if a proof shall be obtained in which all intermediate goals are function-free.) As mentioned, query heads are guaranteed to remain function-free by the correctness of $\mathcal{P}$. For instance, the dependency graph of the goal $\leftarrow a(x)^{(0)},\ b(x,y)^{(1)},\ c(y,z)^{(2)},\ d(z,w)^{(3)}.$ with

$$c_0:\ a(x) \leftarrow a'(x). \qquad c_1:\ b(f(x),x) \leftarrow b'(x).$$
$$c_2:\ c(x,x) \leftarrow c'(x). \qquad c_3:\ d(g(x),x) \leftarrow d'(x).$$

would be $G = \langle \{0,1,2,3\}, \{\langle 1,0\rangle, \langle 3,2\rangle\}\rangle$, i.e. the first subgoal must be resolved before the second and the third subgoal must be resolved before the fourth.

We can now show that such a dependency graph G *is always acyclic*. In fact, if it were not, $\mathcal{P}$ could not be a valid proof, because unification would fail when trying to unify a variable in such a cycle with a function term that *contains that variable*. This is easy to see because each function term given our construction used for obtaining Horn clauses from cind's contains *all* variables appearing in that same (head) atom. Consider for instance

[5] The correctness of the proof $\mathcal{P}$ alone assures that the query head will be function-free as well.

$$q(x) \leftarrow a(x,y),\ a(y,z),\ b(w,z),\ b(z,y).$$

$$\{\langle x,y\rangle \mid \exists z : a(x,z) \wedge a(z,y)\} \supseteq \{\langle x,y\rangle \mid s(x,y)\}$$
$$\{\langle x,y\rangle \mid \exists z : b(x,z) \wedge b(z,y)\} \supseteq \{\langle x,y\rangle \mid s(x,y)\}$$

where s is a source. There is no rewriting under our two semantics, because the dependency graph of our above construction is cyclic already for our initial goal, the body of q.

However, since G is acyclic given a proof $\mathcal{P}$, we can unfold a nonempty set of atoms (those unreachable from other subgoals in graph G) with our intermediate goals until the proof has been completed. □

As an immediate consequence of Lemma 4.2 (which assures that for each resolution proof $\mathcal{P}$ showing $\Sigma \vDash Q \supseteq Q'$ we can produce an equivalent function-free proof $\mathcal{P}'$ that will be covered by Algorithm 4.1), we have

Lemma 4.3. (Completeness of Algorithm 4.1) If $\Sigma \vDash Q \supseteq Q'$ and $Preds(Q') \subseteq \mathbf{S}$, then Algorithm 4.1 computes a conjunctive query Q'' s.t. $\Sigma \vDash Q \supseteq Q'' \supseteq Q'$. □

Lemma 4.1 and Lemma 4.3 taken together imply Theorem 4.2. Let us visualize the implications of Lemma 4.2 with an example.

Example 4.2. Given a boolean conjunctive query $q \leftarrow b(x,x,0).$ and the following set of Horn clauses which, as is easy to see, are the normal form of a set of cind's, which we do not show in order to reduce redundancy.

$b(x',y',0) \leftarrow a(x,y,2),\ e_\epsilon(x,x'),\ e_1(y,y').$ c_0
$b(x',y',2) \leftarrow a(x,y,0),\ e_1(x,x'),\ e_0(y,y').$ c_4, c_{10}, c_{11}
$b(x',y',0) \leftarrow a(x,y,1),\ e_0(x,x'),\ e_\epsilon(y,y').$ c_{12}, c_{18}, c_{19}
$b(x',y',1) \leftarrow a(x,y,0),\ e_1(x,x'),\ e_1(y,y').$ c_{20}, c_{25}

$e_\epsilon(x,x) \leftarrow v(x).$ c_2, c_{17}
$e_1(x,f_1(x)) \leftarrow v(x).$ c_3, c_8, c_{23}, c_{24}
$e_0(x,f_0(x)) \leftarrow v(x).$ c_2, c_{17}
$v(x) \leftarrow b(x,y,s).$ c_5, c_{13}, c_{21}
$v(y) \leftarrow b(x,y,s).$ c_6, c_{14}
$a(x,y,s) \leftarrow b(x,y,s).$ c_1, c_7, c_{15}

where x, y, x', y', s are variables. Let $\mathcal{P}$ be the resolution proof

(0) $\leftarrow b(x,x,0)^{(0)}.$
(1) $\leftarrow a(x,y,2)^{(1)},\ e_\epsilon(x,z)^{(2)},\ e_1(y,z)^{(3)}.$
(2) $\leftarrow b(f_1(y),y,2)^{(4)},\ v(f_1(y))^{(5)},\ v(y)^{(6)}.$
(3) $\leftarrow a(x_1,y_1,0)^{(7)},\ e_1(x_1,f_1(y))^{(8)},\ e_0(y_1,y)^{(9)},\ b(f_1(y),v_1,2)^{(10)},$
$b(v_2,y,2)^{(11)}.$ $\dagger_{10}, \dagger_{11}$
(4) $\leftarrow b(f_0(y_1),y_1,0)^{(12)},\ v(f_0(y_1))^{(13)},\ v(y_1)^{(14)}.$
(5) $\leftarrow a(x_2,y_2,1)^{(15)},\ e_0(x_2,f_0(y_1))^{(16)},\ e_\epsilon(y_2,y_1)^{(17)}, b(f_0(y_1),v_1,0)^{(18)},$
$b(v_2,y_1,0)^{(19)}.$ $\dagger_{18}, \dagger_{19}$
(6) $\leftarrow b(y_1,y_1,1)^{(20)},\ v(y_1)^{(21)}.$
(7) $\leftarrow a(x,x,0)^{(22)},\ e_1(x,f_1(x))^{(23)},\ e_1(x,f_1(x))^{(24)},\ b(y_1,v_1,1)^{(25)}.$ $\dagger_{25}$
(8) $\leftarrow a(x,x,0)^{(22)},\ v(x)^{(26)}.$

which rewrites our query into $q \leftarrow a(x,x,0), v(x).$ and in which we have superscribed each subgoal with its assigned index. In each resolution step, a goal $\leftarrow A^{(i_1)}, \ldots, A^{(i_n)}.$ is unfolded with the clauses $c_{i_1}, \ldots, c_{i_n}$, as annotated above. To keep things short, we have eliminated subgoals (marked with a dagger † and their index) that are redundant with a different branch of the proof. As claimed in our theorem, $\mathcal{P}$ can be transformed into the following proof in which each intermediate step is function-free.

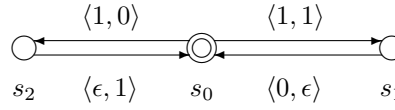
(0) $\leftarrow b(x,x,0)^{(0)}.$
(1) $\leftarrow a(x,y,2)^{(1)},\ e_\epsilon(x,z)^{(2)},\ [e_1(y,z)^{(3)}].$
(2) $\leftarrow b(x,y,2)^{(4)},\ v(x)^{(5)},\ [e_1(y,x)^{(3)}].$
(3) $\leftarrow a(x_1,y_1,0)^{(7)},\ e_1(x_1,x)^{(8)},\ e_0(y_1,y)^{(9)},\ b(x,v_1,2)^{(10)},\ [e_1(y,x)^{(3)}].$ $\dagger_{10}$
(4) $\leftarrow a(x_1,y_1,0)^{(7)},\ e_1(x_1,x)^{(8)},\ [e_0(y_1,y)^{(9)}],\ \not[e_1(y,x)^{(3)}\not].$
(5) $\leftarrow b(y,y_1,0)^{(12)},\ v(y)^{(14)},\ [e_0(y_1,y)^{(9)}].$
(6) $\leftarrow a(x_2,y_2,1)^{(15)},\ e_0(x_2,y)^{(16)},\ e_\epsilon(y_2,y_1)^{(17)},\ b(y,v_1,0)^{(18)},\ \not[e_0(y_1,y)^{(9)}\not].$ $\dagger_{18}$
(7) $\leftarrow b(y_1,y_1,1)^{(20)},\ v(y_1)^{(21)}.$
(8) $\leftarrow a(x_3,y_3,0)^{(22)},\ e_1(x_3,y_1)^{(23)},\ e_1(y_3,y_1)^{(24)},\ b(y_1,v_1,1)^{(25)}.$ $\dagger_{25}$
(9) $\leftarrow a(x_3,x_3,0)^{(22)},\ v(x_3)^{(26)}.$

The subgoals that we have marked with brackets [ ] had been blocked at a certain step to keep the proof function-free. □

Note that Example 4.2 constitutes another encoding of PCP that shows the undecidability of query rewriting with cind's. The PCP instance

$$\mathcal{I} = \{\langle x_1 = 10, y_1 = 1\rangle, \langle x_2 = 1, y_2 = 01\rangle\}$$

is encoded in the first four Horn clauses, which can be viewed as realizing a nondeterministic automaton that accepts two words $x_{i_1} \ldots x_{i_k}$ and $y_{i_1} \ldots y_{i_k}$ if they can be constructed using the dominos of $\mathcal{I}$. In the start state s_0, a domino $\langle x_i, y_i\rangle$ out of $\mathcal{I}$ is chosen. The symbols in x_i and y_i are then accepted one by one. If one of the two words x_i, y_i is longer than the other one, the shorter one is appended ϵ symbols. We return to the state s_0 no sooner than all symbols of a domino have been accepted. For the instance of Example 4.2, we thus have an automaton with three states.



The encoding again allows to show the undecidability of our query rewriting problem (A PCP instance is satisfiable iff the maximally contained rewriting of $q \leftarrow b(x,x,0).$ under Σ is nonempty.) as well as the undecidability of query containment under a set of cind's. (A PCP instance is satisfiable if and only if $\Sigma \vDash \{\langle\rangle \mid \exists x : v(x) \wedge a(x,x,0)\} \subseteq \{\langle\rangle \mid \exists x : b(x,x,0)\}$.)

Of course this correspondence between function-free and general resolution proofs does not hold for Horn clauses in general.

Example 4.3. The boolean query $q \leftarrow a_1(u,v), b_1(u,v).$ and the Horn clauses

$$a_1(f(x), y) \leftarrow a_2(x, y). \qquad a_2(x, g(y)) \leftarrow a_3(x, y).$$
$$b_1(x, g(y)) \leftarrow b_2(x, y). \qquad b_2(f(x), y) \leftarrow b_3(x, y).$$

taken together entail $q \leftarrow a_3(x, y), b_3(x, y).$ even though one cannot arrive at a function-free intermediate rewriting by either unfolding the left subgoal (resulting in $q \leftarrow a_2(x, y),\ b_1(f(x), y).$) or the right subgoal (which would result in $q \leftarrow a_1(x, g(y)),\ b_2(x, y).$) of our query first, neither by unfolding both at once (resulting in $q \leftarrow a_2(x, g(y)),\ b_2(f(x), y).$). □

5 Implementation

Our implementation is based on Algorithm 4.1, but makes use of several optimizations. Directly after parsing, Horn clauses whose head predicates are unreachable from the predicates of the query are filtered out. The same is done with clauses not in the set X computed by

$X := \emptyset$;
do $X := X \cup$
$\{c \in C \mid Preds(Body(c)) \subseteq (Sources \cup \{Pred(Head(c')) \mid c' \in X\})\}$;
while X changed;

We have implemented the simple optimizations known from the Bucket Algorithm [17] and the Inverse Rules Algorithm [11] for answering queries using views which are used to reduce the branching factor in the search process. Beyond that, MiniCon descriptions are computed with an intelligent backtracking method that always chooses to cover subgoals first for which this can be done deterministically (i.e., the number of Horn clauses that are candidates for unfolding with a particular subgoal can be reduced to one), thereby reducing the amount of branching.

In the implementation of the deterministic component of our algorithm for generating MiniCon descriptions, we first check whether the corresponding pairs of terms of two atoms to match unify independently before doing full unification. This allows to detect most violations with very low overhead. Given an appropriate implementation, it is possible to check this property in logarithmic or even constant time.

Our unification algorithm allows to pre-specify variables that may in no case be unified with a function term (e.g., for head variables of queries or atoms already over source predicates). This allows to detect the impossibility to create a function-free rewriting as early as possible.

Every time an MCD m is unfolded with a query to produce an *intermediate* rewriting Q, we compute a query Q' (a partial rewriting) as follows.

$$Body(Q') := \{Body_i(Q) \mid m_i \neq \epsilon\}$$
$$Head(Q') := \langle x_1, \ldots, x_n \rangle \text{ s.t. each } x_i \in Vars(Head(Q)) \cap Vars(Body(Q'))$$

Q' is thus created from the new subgoals of the query that have been introduced using the MCD. If Q' contains non-source predicates, the following check

is performed. We check if our rewriting algorithm produces a nonempty rewriting on Q'. This is carried out in depth-first fashion. If the set of cind's is cyclic, we use a maximum lookahead distance to assure that the search terminates. If Q' is not further rewritable, Q does not need to be further processed but can be dropped. Subsequently, (intermediate) rewritings produced by unfolding queries with MiniCon descriptions are simplified using tableau minimization.

An important performance issue in Algorithm 4.1 is the fact that MCDs are only applied one at a time, which leads to redundant rewritings as e.g. the same MCDs may be applicable in different orders (as is true for the classical problem of answering queries using views, a special case) and thus a search space that may be larger than necessary. We use dependency graph-based optimizations to check if a denser packing of MCDs is possible. For the experiments with layered sets of cind's reported on in Section 6 (Figures 2 and 3), MCDs are packed exactly as densely as in the MiniCon algorithm of [18].

Distribution. The implementation of our query rewriter consists of about 9000 lines of C++ code. Binaries for several platforms as well as examples and a Web demonstrator that allows to run limited-size problems online are available on the Web at [13].

6 Experiments

A number of experiments have been carried out to evaluate the scalability of our implementation. These were executed on a 600 MHz dual Pentium III machine running Linux. A benchmark generator was implemented that randomly generated example chain queries and sets of chain cind's[6]. Chain queries are conjunctive queries of the form

$$q(x_1, x_{n+1}) \leftarrow p_1(x_1, x_2),\ p_2(x_2, x_3),\ \ldots,\ p_{n-1}(x_{n-1}, x_n),\ p_n(x_n, x_{n+1}).$$

Thus, chain queries are constructed by connecting binary predicates via variables to form chains, as shown above. In our experiments, the distinguished (head) variables were the first and the last. The chain cind's had between 3 and 6 subgoals in both the subsuming and the subsumed queries.

In all experiments, the queries had 10 subgoals, and we averaged timings over 50 runs. Sets of cind's were always acyclic. This was ascertained by the use of predicate indexes such that the predicates in a subsumer query of a cind only used indexes greater than or equal to a random number determined for each cind, and subsumed queries only used indexes smaller than that number. Times for parsing the input were excluded from the diagrams, and redundant rewritings were not eliminated[7]. Diagrams relate reasoning times on the (logarithmic-scale) vertical

[6] Experiments with various kinds of random queries and constraints were carried out, too. In this paper, we only report on chain queries, but the experiments with random queries were similarly favorable.

[7] Note that our implementation optionally can make finite rewritings non-redundant and minimal. However, for our experiments, these options were not active.

axis to the problem size expressed by the number of cind's on the horizontal axis.

We report on three experiments.

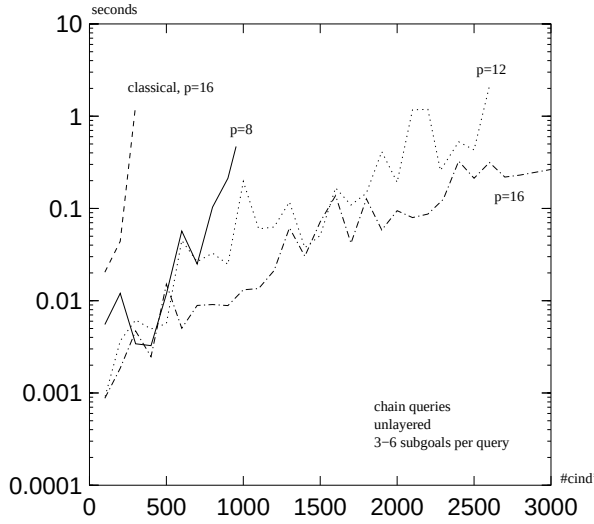


Fig. 1. Experiments with chain queries and nonlayered chain cind's.

Figure 1 shows timings for non-layered sets of constraints. By the steep line on the left we report on an alternative query rewriting algorithm that we have implemented and which follows a traditional resolution strategy. This algorithm (evaluated using instances with 16 predicates) is compared to and clearly outperformed by our new algorithm (with three different numbers of predicates; 8, 12, and 16). Clearly, the more predicates are available, the sparser the constraints get. Thus, more predicates render the query rewriting process simpler.

In Figure 2, we report on the execution times of our new algorithm with cind's generated with an implicit layering[8] of predicates (with 2 layers). This experiment is in principle very similar to local-as-view rewriting with $p/2$ global predicates and $p/2$ source predicates (where the subsumer queries of cind's correspond to logical views in the problem of answering queries using views), followed by view unfolding to account for the subsumed sides of cind's. We again report timings for three different total numbers of predicates, 8, 12, and 16.

In Figure 3, the new algorithm computes maximally contained rewritings for 20 and 40 predicates, which are grouped into stacks of five layers of 4 and 8 predicates each, respectively. Of the five sets of predicates, one constitutes the sources and one the "integration schema" over which queries are asked, and four equally sized sets of cind's bridge between these layers.

[8] See Section 2 for our definition of layered sets of cind's.

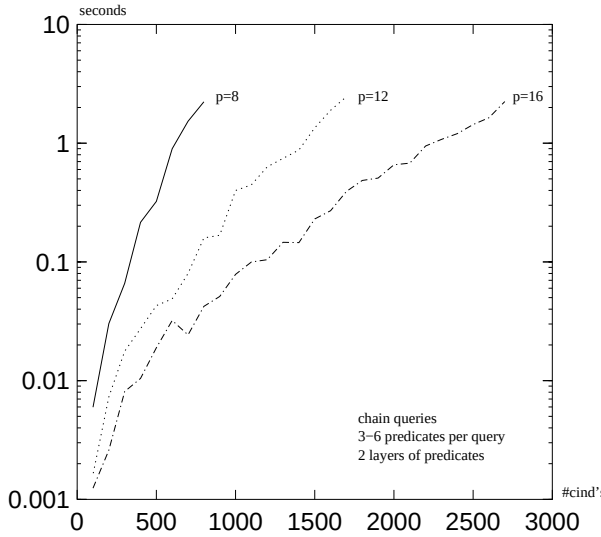


Fig. 2. Experiments with chain queries and two layers of chain cind's.

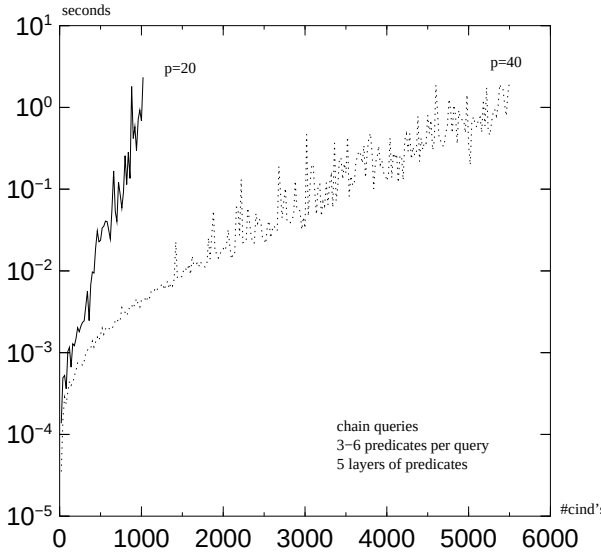


Fig. 3. Experiments with chain queries and five layers of chain cind's.

As can be seen by comparing the second and third diagrams with the first, the hardness of the layered problems is more homogeneous. Particularly in Figure 1 and Figure 2, one can also observe subexponential performance. Note that in the experiment of Figure 3, timings were taken in steps of 20 cind's, while in the other experiments, this step length was 100.

7 Discussion and Conclusions

This paper has addressed the query rewriting problem in data integration from a fresh perspective. We compute maximally contained rewritings with expressive symmetric constraints, which we call Conjunctive Inclusion Dependencies. We have proposed a new query rewriting algorithm based on techniques developed for the problem of answering queries using views (i.e., the MiniCon algorithm), which allows to apply time-tested (e.g., tableau minimization) techniques and algorithms from the database field to the query rewriting problem.

The main advantage of the new algorithm is that intermediate results are (function-free) queries and can be immediately made subject to query optimization techniques. As a consequence, further query rewriting may start from simpler queries, leading to an increase in performance and fewer redundant results that have to be found and later be eliminated. Thus, it is often possible to detect dead ends early. As a trade-off (as can be seen in Algorithm 4.1), an additional degree of nondeterminism is introduced compared to resolution-based algorithms that may temporarily introduce function terms.

In the context of data integration, there are usually a number of regularities in the way constraints are implemented and queries are posed. We expect to have a number of schemata, each one containing a number of predicates. Between the predicates of one schema, no constraints for data integration uses are defined. Moreover, we expect inter-schema constraints to be of the form $Q_1 \subseteq Q_2$ where most (or all) predicates in Q_1 belong to one and the same schema, while the predicates of Q_2 belong to another one. Queries issued against the system are usually formulated in terms of a single schema, and such a layering often propagates along intermediate rewritings. Given these assumptions, we suspect our approach – when optimization techniques from the database area are applied to intermediate results – to have a performance advantage over classical resolution-based algorithms, which do not exploit such techniques.

Our experiments show that our approach scales to very large and complex data integration settings with many schemata.

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